

# Verification of Timed Asynchronous Programs

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## Abstract

In this paper, we address the verification problem for timed asynchronous programs. We associate to each task, a deadline for its execution. We first show that the control state reachability problem for such class of systems is decidable while the configuration reachability problem is undecidable. Then, we consider the subclass of timed asynchronous programs where tasks are always being executed from the same state. For this subclass, we show that the control state reachability problem is PSPACE-complete. Furthermore, we show the decidability for the configuration reachability problem for the subclass.

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## 1 Introduction

One of the well-known design paradigms in concurrent programs is to break a problem into smaller subproblems which are solved asynchronously and concurrently. Each process or thread in the program can then dispatch tasks to other processes, expecting them to be completed by a certain deadline. Each process has a potentially unbounded bag where its pending tasks are stored. In the asynchronous paradigm, one need not wait for time-consuming tasks to be completed to proceed; asynchronous procedure calls are stored in a task buffer, which are executed later, rather than right away. The tasks which are posted asynchronously have deadlines attached to them, and the process or thread, in whose bag the task has been posted, must execute the task within the deadline. In addition to asynchronous procedure calls, one can also make use of synchronous procedure calls where the caller of the procedure blocks until the callee returns. To summarize, an asynchronous program is one that contains procedure calls which are not immediately executed from the calling site, but stored and dispatched in a non-deterministic order by some scheduler(s) at a later point.



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As an example for timed asynchronous programs, we look at *SwingWorker*, an abstract class developed for the Swing library of Java, and is used to perform lengthy GUI interaction tasks in a background thread. While developing applications, sometimes the GUI hangs when it is trying to do some lengthy task. For such purposes, the *SwingWorker* class schedules the execution of this lengthy task on a different thread while the GUI still remains responsive. There are deadlines associated with the background tasks, and if the worker thread which is handling the background task does not finish by the given deadline, then an interrupt is created. To update the user (and GUI) regarding the progress of background tasks, inter-thread communication is allowed.

Writing correct asynchronous programs and reasoning about their correctness is very difficult, since the creation and execution of tasks within deadlines leads to unpredictable behaviours. The verification of asynchronous programs is hence a very challenging topic. A formal model of mutiset pushdown systems for asynchronous recursive programs was presented in [16]. This model consists of a pushdown automaton equipped with a multiset or bag. The automaton adds pending asynchronous method calls to the bag, and the stack executes synchronous recursive method calls. A task can be taken from the bag for execution when the stack is empty. The control state reachability problem was shown to be decidable with an EXPSPACE lower bound under this model. This shows that the case of single-thread asynchronous programs, the reachability problem is very difficult. Subsequently, [8] showed that control state reachability for single-thread asynchronous recursive programs is EXPSPACE-complete. In all these models, time constraints do not play a role in the execution of the asynchronous methods. In the timed setting, [7] considers asynchronous calls of the form `future(p, t)` posted to the task buffer, where  $p$  is a handler and  $t \in \mathbb{N}$ . The idea is that the handler  $p$  will execute the task in  $t$  time units from now. The execution of the program is controlled by logical ticks of a clock. The model proposed in [7] is a generalization of the models in [16] and [11]. [7] shows that safety checking for such programs is undecidable.

The goal of this paper is to investigate the decidability and complexity of the reachability problem for asynchronous non-recursive programs under dense time. We propose a formalism called multiset timed automata (MTA) where each process is modeled as a timed automaton [2]. Each timed automaton is equipped with a bag or multiset. To handle asynchronous method calls, each timed automaton can post a task to the bag of another automaton. These tasks have deadlines attached to them. The deadline is either a natural number  $d \in \mathbb{N}$  or  $\infty$ . When a task is posted to a bag, its age is considered to be 0, and with elapse of time, the age also grows. A task can be executed by the process in whose bag it lies, before the age of the task exceeds the deadline; tasks whose ages have exceeded the deadline will be forever pending. While a main process picks up pending tasks depending on their ages in [7], in our model, a process can execute a pending task in its bag at its will. There are 2 sources of infinity in our model: one coming from dense-time, and the second coming from the unbounded size of the bags of each process. We investigate control state reachability as well as configuration reachability of this model, and show that control state reachability is decidable and EXPSPACE-hard, while configuration reachability is undecidable. We then identify a practically relevant class of MTA where the task execution happens from the same state in each process, and give a PSPACE-complete decision procedure for control state reachability. The configuration reachability also turns out to be decidable for this class.

## Related Work

Most of the existing work (e.g., [3, 4, 6, 8, 11, 13, 14, 16]) on the formal verification of asynchronous programs considers the untimed version. In [7], the authors consider timed constraints on tasks; however, this model is different from the formal model studied in this paper. In

fact, in [7], the authors assume that a task should always be executed by its deadline and the execution of each task is done in logical zero time. In our model, a task whose age has exceeded the deadline will be forever pending. Furthermore, the control state reachability for the model presented in [7] is undecidable while it is decidable for our model. In [5], the authors consider a model similar to the one considered in this paper and show that the coverability problem is decidable using a different technique than ours.

## 2 Preliminaries

In this section, we introduce some notations and definitions that will be used throughout the paper.

### Notations

We use standard notation  $\mathbb{N}$  for the set of naturals, along with  $\infty$ .  $\mathbb{R}$  represents the set of non-negative real numbers. Let  $\mathcal{X}$  be a finite set of variables called *clocks*, taking values from  $\mathbb{R}$ . A *valuation* on  $\mathcal{X}$  is a function  $\nu : \mathcal{X} \rightarrow \mathbb{R}$ . We assume an arbitrary but fixed ordering on the clocks and write  $x_i$  for the  $i$ -th clock. This allows us to treat a valuation  $\nu$  as a vector  $(\nu(x_1), \nu(x_2), \dots, \nu(x_n))$  in  $\mathbb{R}^{|\mathcal{X}|}$ . For a subset of clocks  $X \subseteq \mathcal{X}$  and valuation  $\nu \in \mathbb{R}^{|\mathcal{X}|}$ , we write  $\nu[X:=0]$  for the valuation where  $\nu[X:=0](x) = 0$  if  $x \in X$ , and  $\nu[X:=0](x) = \nu(x)$  otherwise. For  $t \in \mathbb{R}$ , write  $\nu + t$  for the valuation defined by  $\nu(x) + t$  for all  $x \in \mathcal{X}$ . The valuation  $\mathbf{0} \in \mathbb{R}^{|\mathcal{X}|}$  is a special valuation such that  $\mathbf{0}(x) = 0$  for all  $x \in \mathcal{X}$ . For  $a, b \in \mathbb{N}$  and  $a < b$ , the set  $\mathcal{I}$  of time intervals is defined by  $\mathcal{I} := [a, b] \mid [a, a] \mid (a, b] \mid [a, b) \mid (a, b) \mid [a, \infty) \mid (a, \infty)$ . The set of clock constraints, denoted  $\varphi(\mathcal{X})$ , is the set of Boolean formulae over  $\{x \in I \mid x \in \mathcal{X}, I \in \mathcal{I}\}$ . For a constraint  $g \in \varphi(\mathcal{X})$ , and a valuation  $\nu \in \mathbb{R}^{|\mathcal{X}|}$ , we write  $\nu \models g$  to represent the fact that valuation  $\nu$  satisfies the constraint  $g$ . For example,  $(1.1, 0, 10) \models (x_1 \in (0, 2)) \wedge (x_2 \in [0, 0]) \wedge (x_3 \in (1, \infty))$ .

### Timed Automata

Let  $Act$  denote a finite set called actions. A timed automaton (TA) [2] is a tuple  $\mathcal{A} = (L, L^0, Act, \mathcal{X}, E)$  such that (i)  $L$  is a finite set of locations, (ii)  $\mathcal{X}$  is a finite set of clocks, (iii)  $Act$  is a finite alphabet called an action set, (iv)  $E \subseteq L \times \varphi(\mathcal{X}) \times Act \times 2^{\mathcal{X}} \times L$  is a finite set of transitions, and (v)  $L^0 \subseteq L$  is the set of initial locations. A state  $s$  of a timed automaton is a pair  $s = (\ell, \nu) \in L \times \mathbb{R}^{|\mathcal{X}|}$ . A time elapse transition from  $s = (\ell, \nu)$  to  $s' = (\ell', \nu')$  denoted  $s \xrightarrow{t} s'$  is defined iff  $\ell' = \ell$  and  $\nu' = \nu + t$ . Given  $e = (\ell, g, a, Y, \ell') \in E$ , a discrete transition from  $s$  to  $s'$  on  $e$  is written as  $s \xrightarrow{e} s'$ , such that  $\nu \models g$  and  $\nu' = \nu[Y:=0]$ . A run is a finite sequence  $\rho = s_0 \xrightarrow{t_1} s'_0 \xrightarrow{e_1} s_1 \xrightarrow{t_2} s'_1 \xrightarrow{e_2} s_2 \dots s_{n-1} \xrightarrow{t_n} s'_{n-1} \xrightarrow{e_n} s_n$  of states with alternating time elapse transitions and discrete transitions.

### Multisets or Bags

A *multiset* or *bag* over an alphabet  $\Sigma$  is a mapping  $M : \Sigma \mapsto \mathbb{N}$ . For an element  $a \in \Sigma$ , we use  $a \in M$  to denote that  $M(a) \geq 1$ . We use  $\emptyset$  to denote the empty multiset. Given two multisets  $M_1, M_2$  over  $\Sigma$ , we write  $M_1 \leq M_2$  iff  $M_1(a) \leq M_2(a)$  for all  $a \in \Sigma$ .  $M_1 + M_2$  denotes the multiset  $M$  such that  $M(a) = M_1(a) + M_2(a)$  for all  $a \in \Sigma$ . Likewise,  $M_1 - M_2$

denotes, when it is defined (i.e.,  $M_1 \geq M_2$ ), the multiset  $M$  such that  $M(a) = M_1(a) - M_2(a)$  for all  $a \in \Sigma$ . The notation  $M_1 + a$  denotes a multiset  $M_2$  such that  $M_2(a) = M_1(a) + 1$  and  $M_2(b) = M_1(b)$  for all  $b \neq a$ . Likewise,  $M_1 - a$  denotes, when it is defined (i.e.  $M_1(a) \geq 1$ ), a multiset  $M_2$  such that  $M_2(a) = M_1(a) - 1$  and  $M_2(b) = M_1(b)$  for all  $b \neq a$ . The terms multiset and bag will be used interchangeably.

### Timed Petri Nets

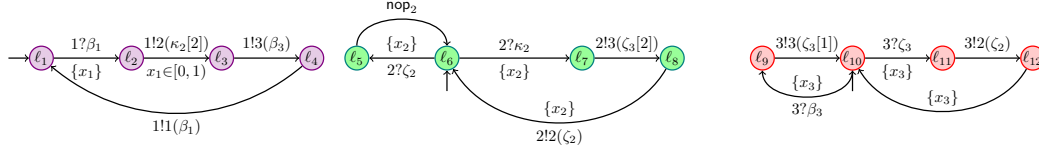
A Timed Petri Net (TPN) [17] is a tuple  $\mathcal{N} = (P, T, F, c)$  where  $P$  is a finite set of places,  $T$  is a finite set of transitions,  $T \cap P = \emptyset$  and  $F \subseteq (P \times T) \cup (T \times P)$  is a flow relation,  $c : F \cap (P \times T) \rightarrow \mathcal{I}$  is a time constraint relation assigning a time interval to every arc from a place to a transition. A marking  $M$  of  $\mathcal{N}$  is a mapping that associates to each place  $p$  a multiset over  $\mathbb{R}$ . A marked TPN is a pair  $(\mathcal{N}, M_0)$  where  $M_0$  is an initial marking, which assigns to each place in  $P$ , an initial multiset of tokens annotated with 0 (the initial age). The dynamics of a TPN consists of two types of transitions rules: firing of a transition and time elapsing. Given  $\mathcal{N}$ , along with a marking  $M$ , (denoted  $(\mathcal{N}, M)$ ) a transition  $t$  is enabled at  $M$  iff for all places  $p$  such that  $(p, t) \in F$ , there exists some  $x \in M(p)$ , and  $x \in c(p, t)$ . If  $t$  is enabled by  $M$ , then it can be fired, producing a marking  $M'$  obtained from  $M$  by (i) removing a token from  $M(p)$  for all places  $p$  such that  $(p, t) \in F$  and whose age satisfies  $c(p, t)$ , and (ii) adding a token with age 0 to  $M(q)$  for all places  $q$  such that  $(t, q) \in F$ . In a time elapse transition, with an elapsing time  $r \in \mathbb{R}$ , the age of all tokens increases by  $r$ . A marked TPN  $(\mathcal{N}, M_0)$  induces a transition system with states are the markings of  $\mathcal{N}$ , and the transition relation consists of time elapsing and firing transitions.

A read arc in a TPN facilitates firing a transition without removing the token. We use  $F^* \subseteq P \times T$  to denote the set of read arcs and  $c^* : F^* \rightarrow \mathcal{I}$  to denote a function that assigns a time interval to each read arc. A transition  $t$  is enabled iff for all places  $p$  such that  $(p, t)^* \in F^*$ , there exists some  $x \in M(p)$  and  $x \in c^*(p, t)$ . The transition system induced by a marked TPN with read arcs can be defined in a similar manner as for marked TPN. A 1-safe marking is one where  $|M(p)| \leq 1$  for all  $p \in P$ . A 1-safe TPN is a marked TPN  $(\mathcal{N}, M_0)$ , with  $F \cap F^* = \emptyset$ , where all markings which are reachable from  $M_0$  are 1-safe.

**Coverability problem.** For markings  $M_1$  and  $M_2$  in a TPN  $\mathcal{N}$ , define  $M_1 \leq M_2$  iff for all  $p \in P$ ,  $M_1(p) \leq M_2(p)$ . The coverability problem for  $\mathcal{N}$  asks whether, given a marking  $M$ , it is possible to reach a marking  $M'$  in  $\mathcal{N}$  from the initial marking  $M_0$  such that  $M \leq M'$ .

## 3 Multiset Timed Automata

Let  $\mathcal{T} = \{T_1, \dots, T_N\}$  be a set consisting of  $N \geq 1$  timed automata  $T_i = (L_i, L_i^0, Act_i, \mathcal{X}_i, E_i)$ . A Multiset Timed Automata (MTA) is defined as  $\mathcal{M} = (\Sigma, \mathcal{T}, \mathcal{X}, St)$ , where  $\Sigma$  is a finite alphabet called *tasks*,  $\mathcal{X} = \biguplus_{i=1}^N \mathcal{X}_i$  is the finite disjoint union of clocks in  $T_i$ ,  $St$  is a function that assigns a finite multiset  $St(i)$  over  $\Sigma$  (possibly empty) to the timed automaton  $T_i$ . This is the initial set of tasks assigned to  $T_i$ . The actions  $Act_i$  are defined as  $Act_i = \{i!j(a[d]), i?a \mid a \in \Sigma, j \in \{1, \dots, N\}, d \in \mathbb{N} \cup \{\infty\}\} \cup \{\text{nop}_i\}$ . The number  $d$  is the *deadline* for the task  $a$ . The action  $i!j(a[d])$  represents  $T_i$  adding the task  $a$  to the bag of automaton  $T_j$ , and the task  $a$  has an associated deadline  $d$ . Likewise, the action  $i?a$  represents automaton  $T_i$  picking up the task  $a$  from its bag, provided its age has not exceeded the deadline. For readability reasons, we assume that any outgoing transition from any initial location is labeled by an action of the form  $i?a$ . We use the notation  $N$ -MTA whenever we need to clarify the number of timed automata  $T_i$  which are used in the definition.



■ **Figure 1** A stateless and time independent 3-MTA consisting of timed automata  $T_1, T_2, T_3$  from left to right. When the deadline of a task is  $\infty$ , we do not mention it.

Let  $\bar{q} = (q_1, \dots, q_N)$  be a tuple of states, where  $q_i = (s_i, \nu_i)$  is the current state of  $T_i$ . Let  $\bar{m} = (M_1, \dots, M_N)$  be a tuple of multisets. Each element in  $M_i$  has the form  $\alpha = (a, r, d) \in \Sigma \times \mathbb{R} \times \mathbb{N}$  consisting of *pending tasks*, their *ages*, and their *deadlines* in  $T_i$ . The age of a task in a bag is the time elapse since it has been added to the bag. For  $t \in \mathbb{R}$ , let  $\bar{q} + t$  represent the tuple  $(q'_1, \dots, q'_N)$  where  $q'_i = (s_i, \nu_i + t)$ . For an element  $\alpha = (a, r, d) \in M_i$ ,  $\alpha + t = (a, r + t, d)$ ;  $M_i + t$  is the multiset obtained by replacing each  $\alpha \in M_i$  with  $\alpha + t$ . We define  $\bar{m} + t$  as the tuple  $(M_1 + t, \dots, M_N + t)$ .

A configuration  $\mathbf{c}$  of an  $N$ -MTA is the tuple  $(\bar{q}, \bar{m})$  consisting of the current states of all the  $N$  timed automata, along with the multisets of pending tasks corresponding to each  $T_i$ . An initial configuration is defined as  $\mathbf{c}_0 = (\bar{q}_0, \bar{m}_0)$ , where  $\bar{q}_0$  is the tuple  $((\ell_1^0, \mathbf{0}), \dots, (\ell_N^0, \mathbf{0}))$  of initial states of all  $T_i$  ( $\ell_i^0 \in L_i^0$ ) and  $\bar{m}_0 = (M_1, \dots, M_N)$  where  $M_i((a, r, d)) = St(i)(a)$ , for all  $a \in \Sigma$ ,  $r = 0$  and  $d = \infty$ , and  $M_i((a, r, d)) = 0$  otherwise. Given two configurations  $\mathbf{c} = (\bar{q}, \bar{m})$ , and  $\mathbf{c}' = (\bar{q}', \bar{m}')$ , we have:

- For  $t \in \mathbb{R}$ ,  $\mathbf{c} \xrightarrow{t} \mathbf{c}'$  is a time elapse transition iff  $\bar{q}' = \bar{q} + t$ ,  $\bar{m}' = \bar{m} + t$ .
- Let  $e_i = (\ell_i, g_i, act_i, Y_i, \ell'_i) \in E_i$ . Then,  $\mathbf{c} \xrightarrow{e_i} \mathbf{c}'$  iff
  - $q_i = (\ell_i, \nu_i)$ ,  $\nu_i \models g_i$ ,  $q'_i = (\ell'_i, \nu'_i)$ ,  $\nu'_i = \nu_i[Y_i := 0]$ , and for all  $k \neq i$ ,  $q'_k = q_k$ , and,
  - If  $act_i = i!j(a[d])$ , then  $M'_j = M_j + (a, 0, d)$ , and  $M'_k = M_k$  for all  $k \neq j$ ,
  - If  $act_i = i?a$ , then  $\exists c, d$ , such that  $(a, c, d) \in M_i$ ,  $M'_i = M_i - (a, c, d)$ , and  $c \leq d$  (i.e. the age of the task has not yet exceeded the deadline) and  $M'_k = M_k$  for all  $k \neq i$ ,
  - If  $act_i = \text{nop}_i$ , then  $M'_k = M_k$  for all  $1 \leq k \leq N$ .

Starting with an initial configuration  $\mathbf{c}_0$ , a run  $\rho$  is defined as a finite sequence of alternating time elapse and discrete transitions of the form  $\mathbf{c}_0 \xrightarrow{t_0} \mathbf{c}'_0 \xrightarrow{e_1} \mathbf{c}_1 \xrightarrow{t_1} \mathbf{c}'_1 \xrightarrow{e_2} \mathbf{c}_2 \dots \xrightarrow{e_j} \mathbf{c}_j$  or  $\mathbf{c}_0 \xrightarrow{t_0} \mathbf{c}'_0 \xrightarrow{e_1} \mathbf{c}_1 \xrightarrow{t_1} \mathbf{c}'_1 \xrightarrow{e_2} \mathbf{c}_2 \dots \xrightarrow{t_j} \mathbf{c}'_j$ . In that case we say the configuration  $\mathbf{c}_j$  is reachable from the initial configuration  $\mathbf{c}_0$  by the run  $\rho$ .

In this paper, we consider the following problems. Let  $\bar{s} = (s_1, \dots, s_N) \in L_1 \times \dots \times L_N$ .

- P1 Control State Reachability.** Given a particular tuple of locations  $\bar{s} = (s_1, \dots, s_N)$  of an  $N$ -MTA  $\mathcal{M}$ , the control state reachability problem asks if starting from the initial configuration  $\mathbf{c}_0$  of  $\mathcal{M}$ , there is a run reaching a configuration  $\mathbf{c} = (\bar{q}, \bar{m})$  such that  $q_i = (s_i, \nu_i)$  for some  $\bar{m}$ , and for some  $\nu_i$ , for all  $1 \leq i \leq N$ .
- P2 Configuration Reachability.** Given a particular tuple of locations  $\bar{s} = (s_1, \dots, s_N)$  of an  $N$ -MTA  $\mathcal{M}$ , the configuration reachability problem asks if starting from the initial configuration  $\mathbf{c}_0$  of  $\mathcal{M}$ , there is a run reaching a configuration  $\mathbf{c} = (\bar{q}, \bar{m})$  such that  $\bar{m} = (\emptyset, \dots, \emptyset)$  and  $q_i = (s_i, \mathbf{0})$ , for all  $1 \leq i \leq N$ .

### Stateless and Time-Independent MTA

An  $N$ -MTA is said to be *stateless* if  $E_i \cap (L_i \setminus \{\ell_i^0\} \times \varphi(\mathcal{X}_i) \times \{i?a | a \in \Sigma\} \times 2^{\mathcal{X}_i} \times L_i) = \emptyset$  for all  $1 \leq i \leq N$ , and some  $\ell_i^0 \in L_i^0$ . The stateless condition ensures that a new task can be picked by an automaton only from a unique initial location. An  $N$ -MTA is said to be

*time-independent* iff, in each  $T_i$ , all clocks are reset on picking a task from the multiset, and no clock constraints are checked (i.e.  $E_i \cap (L_i \times \varphi(\mathcal{X}_i) \times \{i?a|a \in \Sigma\} \times (2^{\mathcal{X}_i} \setminus \{\mathcal{X}_i\}) \times L_i) = \emptyset$  and  $E_i \cap (L_i \times (\varphi(\mathcal{X}_i) \setminus \{true\}) \times \{i?a|a \in \Sigma\} \times 2^{\mathcal{X}_i} \times L_i) = \emptyset$  for all  $1 \leq i \leq N$ ).

Figure 1 describes a stateless and time-independent MTA  $\mathcal{M}$  consisting of 3 timed automata  $T_1, T_2, T_3$ . The following is a run in  $\mathcal{M}$ . The initial configuration  $\mathbf{c}_0 = (\bar{q}_0, \bar{m}_0)$  where  $\bar{q}_0 = ((\ell_1, 0), (\ell_6, 0), (\ell_{10}, 0))$  and  $\bar{m}_0 = (M_1, M_2, M_3)$  with multisets  $M_1 = \{(\beta_1, 0, \infty)\}$ ,  $M_2 = \{(\beta_2, 0, \infty)\}$ , and  $M_3 = \{(\beta_3, 0, \infty)\}$ . Let  $e_{i,j}$  denote the transition from location  $\ell_i$  to  $\ell_j$  (in the example, we have at most one transition between any pair of locations  $\ell_i, \ell_j$ ). For example,  $e_{23} = (\ell_2, x_1 \in [0, 1], 1!2(\kappa_2[2]), \emptyset, \ell_3)$ . Consider the run  $\sigma$   $\mathbf{c}_0 \xrightarrow{0.5} \mathbf{c}'_0 \xrightarrow{e_{1,2}} \mathbf{c}_1 \xrightarrow{0.3} \mathbf{c}'_1 \xrightarrow{e_{2,3}} \mathbf{c}_2 \xrightarrow{0.5} \mathbf{c}'_2 \xrightarrow{e_{6,7}} \mathbf{c}_3 \xrightarrow{0.2} \mathbf{c}'_3 \xrightarrow{e_{7,8}} \mathbf{c}_4 \xrightarrow{0} \mathbf{c}'_4 \xrightarrow{e_{10,9}} \mathbf{c}_5 \xrightarrow{0.4} \mathbf{c}'_5 \xrightarrow{e_{3,4}} \mathbf{c}_6 \xrightarrow{0.1} \mathbf{c}'_6 \xrightarrow{e_{4,1}} \mathbf{c}_7 \xrightarrow{0.1} \mathbf{c}'_7 \xrightarrow{e_{1,2}} \mathbf{c}_8 \xrightarrow{0.1} \mathbf{c}'_8 \xrightarrow{e_{8,6}} \mathbf{c}_9 \xrightarrow{0.2} \mathbf{c}'_9 \xrightarrow{e_{9,10}} \mathbf{c}_{10} \xrightarrow{0.6} \mathbf{c}'_{10} \xrightarrow{e_{6,5}} \mathbf{c}_{11} \xrightarrow{0.5} \mathbf{c}'_{11} \xrightarrow{e_{10,11}} \mathbf{c}_{12} \xrightarrow{0.9} \mathbf{c}'_{12} \xrightarrow{e_{11,12}} \mathbf{c}_{13}$  which reaches locations  $(\ell_2, \ell_5, \ell_{12})$  in  $T_1, T_2, T_3$  respectively.

## 4 Control State Reachability

In the following, we first prove that the control reachability is decidable with a non-primitive complexity (at the level  $F_{\omega^\omega}$  in the fast growing hierarchy [9]). Then, we show that the control state reachability for stateless and time independent MTA is PSPACE-complete.

► **Theorem 1.** *The control state reachability problem for  $N$ -MTA is reducible to the coverability problem for timed Petri nets with read-arcs.*

**Proof.** We give a translation from an  $N$ -MTA  $\mathcal{M}$  to a TPN with read arcs  $\mathcal{N}$  such that the control state reachability of  $\mathcal{M}$  reduces to the coverability of  $\mathcal{N}$ .

Let  $\mathcal{M} = (\Sigma, \mathcal{T}, \mathcal{X}, St)$  be an  $N$ -MTA. Without loss of generality, assume that we are interested in reaching  $\bar{f} = (f_1, \dots, f_N) \in L_1 \times \dots \times L_N$ . Given the  $N$ -MTA  $\mathcal{M}$ , we construct a timed Petri net  $\mathcal{N}$  as follows. There is a place  $p_\ell$  in the net corresponding to each location  $\ell \in L_i$  in  $T_i$  for each  $i \in \{1, \dots, N\}$ . For each  $T_i$ , there is one and only one marked place  $p_\ell$  such that  $\ell \in L_i$ , to denote that the control of  $T_i$  is at a certain location  $\ell$ . For each clock  $x$  in  $\mathcal{X}$ , we have a place  $p_x$  in the net. Next, we model the multisets  $M_i$  of each  $T_i$ . Let  $d_{max} \in \mathbb{N}$  be the maximal value used for any deadline in  $\mathcal{M}$ . The possible task, deadline combinations are in the set  $\Sigma \times \{0, 1, \dots, d_{max}, \infty\}$ . Therefore, corresponding to each  $T_i$ , we have  $|\Sigma| \times (d_{max} + 2)$  places in the net. We need to have these many places so as to distinguish between the tokens. Thus, for each pair  $(a, d) \in \Sigma \times \{0, 1, \dots, d_{max}, \infty\}$ , we have the places  $p_{(a,d)}^1, \dots, p_{(a,d)}^N$ .

A transition of the form  $(\ell, g, i?a, Y, \ell')$  in automaton  $T_i$  is simulated by a transition in  $\mathcal{N}$  as follows. A token from the place  $p_\ell$  corresponding to the location  $\ell$ , and a token from one of the places  $p_{(a,d)}^i, d \in \{0, 1, \dots, d_{max}, \infty\}$  are removed. The deadline is checked on the arc via a constraint  $[0, z]$  from the place  $p_{(a,z)}^i$  containing the token. A token is added to the place  $p_{\ell'}$  corresponding to  $\ell'$ . A transition of the form  $(\ell, g, i!j(a[d]), Y, \ell')$  in automaton  $T_i$  is simulated in a similar way. The tokens corresponding to locations are removed and added as in the previous case and a token is added to the place  $p_{(a,d)}^j$ . The clock constraints corresponding to any transition are checked using read arcs from the places simulating the clocks. Clock resets are simulated by removing a token and putting back a token in the place corresponding to the clock.

The details of the formal construction of  $\mathcal{N}$  and the correctness proof can be found in the extended version of the paper [1]. ◀

As a corollary of Theorem 1, we get:

► **Corollary 2.** *The control state reachability problem for (time-independent)  $N$ -MTA is decidable.*

Observe that we can easily show that the coverability of Petri nets is reducible to the control state reachability problem for time-independent (resp. stateless)  $N$ -MTA (in the same way as the proof of EXPSPACE lower bound for the model multiset pushdown systems presented in [16]). Therefore, the control state reachability problem for time-independent (resp. stateless)  $N$ -MTA is EXPSPACE-hard.

In the rest of this section, we consider the case of stateless and time-independent  $N$ -MTA.

► **Theorem 3.** *The control state reachability problem for stateless and time-independent  $N$ -MTA is PSPACE-complete (for  $N \geq 1$ ).*

**Proof.** Since MTA subsume timed automata [2], the PSPACE-hardness of the control state reachability of MTA follows directly from the PSPACE-hardness of reachability of timed automata. The rest of the proof is devoted to proving the PSPACE-membership of the problem.

Let  $\mathcal{M} = (\Sigma, \mathcal{T}, \mathcal{X}, St)$  be a stateless, time-independent  $N$ -MTA, with  $\mathcal{T} = \{T_1, \dots, T_N\}$ ,  $\mathcal{X} = \bigsqcup_{i=1}^N \mathcal{X}_i$  and  $St$ , the function that assigns an initial multiset  $St(i)$  to each timed automaton  $T_i$ . Incurring a polynomial blowup in the size, we give a reduction from the control state reachability of  $\mathcal{M}$  to the coverability in 1-safe timed Petri net with read arcs. The coverability of 1-safe timed Petri nets with read arcs is known to be PSPACE-complete [17] and our result follows from this.

Without loss of generality, assume that we are interested in reaching  $\bar{f} = (f_1, \dots, f_N) \in L_1 \times \dots \times L_N$ . Let  $\sigma$  be any run from the initial configuration  $\mathbf{c}_0$  of  $\mathcal{M}$  which leads into a configuration with locations  $\bar{f}$ . Let  $\mathbf{c} = (\bar{q}, \bar{m})$  be any configuration that appears in  $\sigma$ . Our proof is divided into two parts.

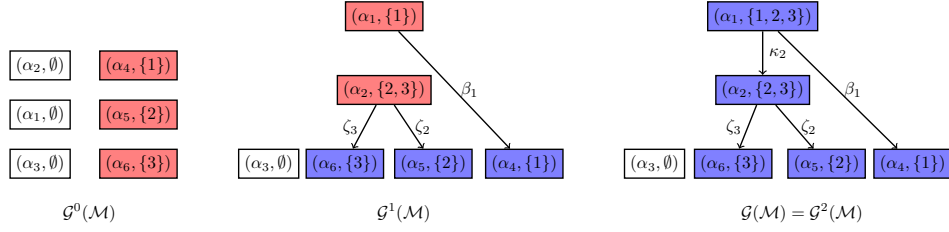
1. We show that the number of *relevant* task tuples along  $\sigma$  is bounded by  $N$ . Intuitively, A task tuple  $(a, r, d) \in \Sigma \times \mathbb{R} \times \mathbb{N}$  is relevant for an automaton  $T_i$  if  $(a, r, d) \in M_j$ , for some  $j$ ,  $(r \leq d)$  and the task  $(a, r, d)$  must be executed by  $T_j$  in order to reach the location  $f_i$ . The irrelevant task tuples can hence be ignored from each  $M_i$ , as they do not affect the control state reachability.
2. The bound on the number of relevant task tuples obtained in the previous step is used in constructing a reachability preserving 1-safe timed Petri net with read arcs.

### Bounding the number of relevant task tuples

Consider the run  $\sigma$  as described above. Starting from  $\mathbf{c}_0$ , let  $\sigma_i$  denote the sequence of transitions (in the order they appear in  $\sigma$ ), pertaining only to  $T_i$ . In the run  $\sigma$  pertaining to the example in figure 1,  $\sigma_1$  consists of all the violet discrete transitions separated by time elapses:  $\mathbf{c}_0 \xrightarrow{0.5} \mathbf{c}'_0 \xrightarrow{e_{1,2}} \mathbf{c}_1 \xrightarrow{0.3} \mathbf{c}'_1 \xrightarrow{e_{2,3}} \mathbf{c}_2 \xrightarrow{1.1} \mathbf{c}'_5 \xrightarrow{e_{3,4}} \mathbf{c}_6 \xrightarrow{0.1} \mathbf{c}'_6 \xrightarrow{e_{4,1}} \mathbf{c}_7 \xrightarrow{0.1} \mathbf{c}'_7 \xrightarrow{e_{1,2}} \mathbf{c}_8$ . We now define a block.

A *block* in  $\sigma_i$  begins with a discrete transition of the form  $\xrightarrow{i?a}$  (for some task  $a$ ) and extends until the next transition of the form  $\xrightarrow{i?b}$  (for some task  $b$ ) is encountered. Thus, a block is a sequence of transitions between two executions of tasks by some  $T_i$ , and begins with some task execution.  $\sigma_1$  has two blocks: the sequence of transitions from  $\mathbf{c}'_0$  till  $\mathbf{c}'_7$  forms a block, and the second block is the transition from  $\mathbf{c}'_7$  to  $\mathbf{c}_8$ . Omitting the time elapse transitions in  $\sigma_i$ , let us label each transition in  $\sigma_i$  with a unique name. Doing this for all  $\sigma_i$  gives us a unique label for each discrete transition in  $\sigma$ . Let  $\mathcal{L} = \{\alpha_1, \dots, \alpha_m\}$  be the set of block labels occurring in  $\sigma$ . In our running example, using labels  $\{\alpha_1, \dots, \alpha_6\}$ , we can label the blocks in  $\sigma$  as  $\mathbf{c}_0 \xrightarrow{0.5} \mathbf{c}'_0 \xrightarrow{e_{1,2}, \alpha_1} \mathbf{c}_1 \xrightarrow{0.3} \mathbf{c}'_1 \xrightarrow{e_{2,3}, \alpha_1} \mathbf{c}_2 \xrightarrow{0.5} \mathbf{c}'_2 \xrightarrow{e_{6,7}, \alpha_2} \mathbf{c}_3 \xrightarrow{0.2} \mathbf{c}'_3 \xrightarrow{e_{7,8}, \alpha_2} \mathbf{c}_4 \xrightarrow{0} \mathbf{c}'_4 \xrightarrow{e_{10,9}, \alpha_3} \mathbf{c}_5 \xrightarrow{0.4} \mathbf{c}'_5 \xrightarrow{e_{3,4}, \alpha_1} \mathbf{c}_6 \xrightarrow{0.1} \mathbf{c}'_6 \xrightarrow{e_{4,1}, \alpha_1} \mathbf{c}_7 \xrightarrow{0.1} \mathbf{c}'_7 \xrightarrow{e_{1,2}, \alpha_4} \mathbf{c}_8 \xrightarrow{0.1} \mathbf{c}'_8 \xrightarrow{e_{8,6}, \alpha_2} \mathbf{c}_9 \xrightarrow{0.2} \mathbf{c}'_9 \xrightarrow{e_{9,10}, \alpha_3} \mathbf{c}_{10} \xrightarrow{0.6} \mathbf{c}'_{10} \xrightarrow{e_{6,5}, \alpha_5} \mathbf{c}_{11} \xrightarrow{0.5} \mathbf{c}'_{11} \xrightarrow{e_{10,11}, \alpha_6} \mathbf{c}_{12} \xrightarrow{0.9} \mathbf{c}'_{12} \xrightarrow{e_{11,12}, \alpha_6} \mathbf{c}_{13}$ . From here on, we refer to the blocks using the block labels.





■ **Figure 2** The dependency graph in stages.  $\mathcal{G}^0(\mathcal{M})$  is the initial graph with no edges.  $\mathcal{G}^{i+1}(\mathcal{M})$  is obtained from  $\mathcal{G}^i(\mathcal{M})$  by changing the color of all the red vertices in  $\mathcal{G}^i(\mathcal{M})$ . The graph stabilizes when there are no red vertices.

For each timed automaton  $T_i$ , we now analyze the blocks which contribute in reaching the desired location  $f_i$ . The last block  $\alpha$  of  $\sigma_i$ , which contains the last task tuple  $(a, r, d)$  executed by  $T_i$  definitely contributes to  $T_i$  reaching  $f_i$ . Likewise, the block  $\alpha'$  which added this last task  $a$  to the bag of  $T_i$  also contributes to  $T_i$  reaching  $f_i$  (note that block  $\alpha'$  may start with a task  $b$  which is executed by  $T_j$ ,  $j \neq i$ ). We can continue backwards in this manner and say that the block  $\alpha''$  which added the task  $b$  to the bag of  $T_j$  also contributes to  $T_i$  reaching  $f_i$  and so on. Given a block label  $\alpha$ , let  $\text{dep}(\alpha)$  denote the set of timed automata  $T_i$  such that  $\alpha$  contributes to  $T_i$  reaching  $f_i$ . Thus, if  $\alpha$  is the last block in  $T_i$ , then  $i \in \text{dep}(\alpha)$  (we just write the indices  $i$  rather than  $T_i$ ). Likewise, if  $i \in \text{dep}(\alpha)$  and if  $\alpha'$  is the block which added the task  $a$  which was executed at the beginning of  $\alpha$ , then  $i \in \text{dep}(\alpha')$ , and so on.  $\text{dep}(\alpha)$  is called the dependency set of  $\alpha$ . In our running example above,  $3 \in \text{dep}(\alpha_6)$  since  $\alpha_6$  is the last block for  $T_3$ ; however the task  $\zeta_3$  which was executed in block  $\alpha_6$  was added in block  $\alpha_2$  ( $e_{7,8}$ ), and the task  $\kappa_2$  executed in block  $\alpha_2$  was added in block  $\alpha_1$ . Thus,  $3 \in \text{dep}(\alpha_6), \text{dep}(\alpha_2), \text{dep}(\alpha_1)$ .

We construct a *dependency graph*  $\mathcal{G}(\mathcal{M})$  which keeps track of the dependencies between blocks. Define a function  $g : \mathcal{L} \rightarrow (\Sigma \times \{1, \dots, N\} \times \mathcal{L}) \cup \{\perp\}$  which maps a block label  $\alpha$  to the triple  $(a, i, \alpha')$  if block  $\alpha$  begins with  $(i?a)$  the execution of task  $a$ , which was added to the bag of  $T_i$  by block  $\alpha'$ . If  $a$  is part of the initial multiset ( $a \in \text{St}(i)$ ) then  $g(\alpha) = \perp$ . The vertex set of  $\mathcal{G}(\mathcal{M})$  is the set of pairs  $(\alpha, \text{dep}(\alpha))$  where  $\alpha$  is a block label and  $\text{dep}(\alpha)$  is its dependency set.  $\mathcal{G}(\mathcal{M})$  is a graph with colored vertices, and is built inductively. To begin, there are no edges, and we have the following vertices.

- Vertices  $(\alpha, \{i\})$  and  $\alpha$  is the last block of  $T_i$ . To begin, we are sure of  $i \in \text{dep}(\alpha)$ . We color these vertices red.
- Vertices  $(\alpha, \emptyset)$ , and  $\alpha$  is not the last block for any  $T_i$ . To begin, we have not yet discovered whether  $\alpha$  contributes to any  $T_i$ , so we keep  $\text{dep}(\alpha) = \emptyset$ . The information with respect to  $\text{dep}(\alpha)$  will be updated when we discover that  $\alpha$  contributed to some  $T_i$ . We color these vertices white.

To add the edges, we repeat the following procedure until no red vertices remain. In each step, we choose a red vertex  $(\alpha, \text{dep}(\alpha))$  and do the following.

1. If  $g(\alpha) = \perp$ , then color  $(\alpha, \text{dep}(\alpha))$  blue,
2. If  $g(\alpha) = (a, i, \alpha')$  and  $(\alpha', \text{dep}(\alpha'))$  is white, then color  $(\alpha, \text{dep}(\alpha))$  blue and color  $(\alpha', \text{dep}(\alpha'))$  red. Update  $\text{dep}(\alpha')$  to be  $\text{dep}(\alpha') \cup \text{dep}(\alpha)$ , and add an edge  $\xrightarrow{a}$  from  $(\alpha', \text{dep}(\alpha'))$  to  $(\alpha, \text{dep}(\alpha))$ .
3. If  $g(\alpha) = (a, i, \alpha')$  and  $(\alpha', \text{dep}(\alpha'))$  is not white, then color  $(\alpha, \text{dep}(\alpha))$  blue, update  $\text{dep}(\alpha')$  to be  $\text{dep}(\alpha') \cup \text{dep}(\alpha)$ , and add an edge  $\xrightarrow{a}$  from  $(\alpha', \text{dep}(\alpha'))$  to  $(\alpha, \text{dep}(\alpha))$ .



Finally, we update the dependency relation  $\text{dep}$  of the vertices as follows: If  $(\alpha, \text{dep}(\alpha))$  and  $(\alpha', \text{dep}(\alpha'))$  are blue with  $g(\alpha) = (a, i, \alpha')$ , then update  $\text{dep}(\alpha')$  to be  $\text{dep}(\alpha') \cup \text{dep}(\alpha)$ . Note that the above procedure terminates, since the number of blue vertices in each step increases. The final graph obtained as result is  $\mathcal{G}(\mathcal{M})$ . Figure 2 describes constructing  $\mathcal{G}(\mathcal{M})$  for the run  $\rho$  discussed above. Consider any vertex  $(\alpha, \text{dep}(\alpha))$  colored blue in  $\mathcal{G}(\mathcal{M})$ . Clearly, this vertex contributes to all  $T_i$  such that  $i \in \text{dep}(\alpha)$ . Consider any path in  $\mathcal{G}(\mathcal{M})$  from a vertex with no incoming edges to a vertex with no outgoing edges. There is at least one such path since the last task executed along  $\sigma$  corresponds to the last block of some  $T_i$  which has not contributed to any  $T_j$ . A path  $v_1 \dots v_s$  in  $\mathcal{G}(\mathcal{M})$  is a dependency path for automaton  $T_i$  if the vertex  $v_s = (\alpha, \text{dep}(\alpha))$ , and  $\alpha$  is the last block for  $T_i$ . Let us go back to our running example run  $\sigma$  using Figure 2. The tasks appearing on the edges of  $\mathcal{G}(\mathcal{M})$  are the relevant tasks. From  $\mathcal{G}(\mathcal{M})$ , the relevant task in the bags when the second block  $\alpha_2$  started is  $\kappa_2$ .  $\kappa_2$  is executed at the beginning of block  $\alpha_2$ . Relevant tasks  $\zeta_2, \zeta_3$  are added to the bag in block  $\alpha_2$ , and  $\alpha_1$  adds  $\beta_1$ .  $\beta_1$  is executed in block  $\alpha_4$  while  $\zeta_2, \zeta_3$  respectively are executed in blocks  $\alpha_5, \alpha_6$ . The relevant tasks along run  $\sigma$  are  $\beta_1, \kappa_2, \zeta_2, \zeta_3$ , of which at most 3 are stored across bags at any point of time. Thus, we can obtain another run  $\sigma'$  which is reachability equivalent to  $\sigma$  as follows. The block  $\alpha_3$  is useless as it is not contributing to any of the automata. Each block begins at a unique initial location of some automaton, and, on the transition which executes the task, it does not check any constraints, and resets all clocks on the transition. Due to this, we can “prune away” a block from a run, and reconnect the run at a later block if we maintain the time elapse in the interim. Hence, removing a useless block of some automaton  $T_i$  does not affect the control reachability, since the next useful block of  $T_i$  again starts from the same initial location of  $T_i$ . Accounting for the time elapse in the useless block is sufficient to ensure that the ages of the pending tasks are accurate.  $\sigma'$  can be constructed with the rest of the blocks, using only the relevant tasks:  $c_0 \xrightarrow{0.5} c'_0 \xrightarrow{e_{1,2}, \alpha_1} c_1 \xrightarrow{0.3} c'_1 \xrightarrow{e_{2,3}, \alpha_1} c_2 \xrightarrow{0.5} c'_2 \xrightarrow{e_{6,7}, \alpha_2} c_3 \xrightarrow{0.2} c'_3 \xrightarrow{e_{7,8}, \alpha_2} c_4 \xrightarrow{0} c'_4 \xrightarrow{0.4} c'_5 \xrightarrow{e_{3,4}, \alpha_1} c_6 \xrightarrow{0.1} c'_6 \xrightarrow{e_{4,1}, \alpha_1} c_7 \xrightarrow{0.1} c'_7 \xrightarrow{e_{1,2}, \alpha_4} c_8 \xrightarrow{0.1} c'_8 \xrightarrow{e_{8,6}, \alpha_2} c_9 \xrightarrow{0.2} c'_9 \xrightarrow{0.6} c_{10} \xrightarrow{e_{6,5}, \alpha_5} c_{11} \xrightarrow{0.5} c'_{11} \xrightarrow{e_{10,11}, \alpha_6} c_{12} \xrightarrow{0.9} c'_{12} \xrightarrow{e_{11,12}, \alpha_6} c_{13}$ .

We want to prove that in any configuration  $\mathbf{c} = (\bar{q}, \bar{m})$  appearing in the run  $\sigma$ , the number of pending tasks maintained in  $\bar{m} = (M_1, \dots, M_N)$  which contribute, in reaching the desired control states, in  $\sigma$  is  $\leq N$ . These are the relevant tasks, and each one is part of a block  $\alpha$ , and the corresponding vertex  $(\alpha, \text{dep}(\alpha))$  in  $\mathcal{G}(\mathcal{M})$  is colored blue. If we attach the color of the vertex  $(\alpha, \text{dep}(\alpha))$  to the task  $a$  in  $g(\alpha)$ , then we want to prove that in any configuration appearing in  $\sigma$ , the number of blue tasks is  $\leq N$ . Assume that there is some configuration  $\mathbf{c} = (\bar{q}, \bar{m})$  in  $\sigma$  such that the number of blue tasks in  $\bar{m}$  is  $p > N$ . Let  $a_1, \dots, a_p$  be the tasks in  $\bar{m}$ , and let  $\alpha_1, \dots, \alpha_p$  be the blocks where these are executed. Since  $p > N$ , and there are only  $N$  multisets in  $\bar{m}$ , there are at least two tasks  $a_i, a_j$  such that  $\text{dep}(\alpha_i) \cap \text{dep}(\alpha_j) \neq \emptyset$ . Observe that, by definition, we have  $\text{dep}(\alpha_k) \neq \emptyset$  for all  $k \in \{1, \dots, p\}$ . Let us assume that  $k \in \text{dep}(\alpha_i) \cap \text{dep}(\alpha_j)$ . Since both are blue, both get executed in  $\sigma$ , and both lie in the dependency path of the last block of the automaton  $T_k$ . Clearly, one must come before the other, and the earlier block has contributed to the creation of the later block. Hence, they cannot be pending at the same time. Thus, the number of blue pending tasks in any configuration is bounded above by  $N$ .

### Construction of 1-safe TPN with read arcs

Now, we are ready to propose a 1-safe timed Petri net (with read-arcs) whose coverability problem is equivalent to the control state reachability problem of the given  $N$ -MTA.

Given the  $N$ -MTA  $\mathcal{M}$  consisting of timed automata  $T_1, \dots, T_N$ , we construct a 1-safe TPN  $\mathcal{N}$ . There is a place  $p_\ell$  corresponding to each location  $\ell \in L_i$  in  $T_i$ . For each  $T_i$ , there is one and only one marked place  $p_\ell$  at any point in the execution, such that  $\ell \in L_i$ , to denote that the control of  $T_i$  is at a certain location  $\ell$ . For each clock  $x$  in  $X$ , there is a place  $p_x$ . Next, we model the multisets  $M_i$  of each  $T_i$ . Let  $d_{max} \in \mathbb{N}$  be the maximal value used for any deadline in  $\mathcal{M}$ . For each task  $a \in \Sigma$ , we have  $|\Sigma| \times (2 + d_{max})$  possible combinations of tasks and associated deadlines. The bound established above tells us that there are at most  $N$  pending tasks in any configuration i.e. at any point we will have to keep track of  $N$  tasks but they can be distributed in any of the multisets. There are  $|\Sigma| \times (d_{max} + 2)$  possibilities for task, deadline pairs. Tasks will be modeled as tokens in the net. So to be able to distinguish between them, for each  $T_i$ , we need  $N \times |\Sigma| \times (d_{max} + 2)$  places ( $N$ , because 1-safe). For each  $T_i$  and for each pair  $(a, d) \in \Sigma \times \{0, 1, \dots, d_{max}, \infty\}$ , we have  $N$  places  $p_{(a,d,1)}^i, \dots, p_{(a,d,N)}^i$ .

A transition of the form  $(\ell_i^0, g, i?a, Y, \ell')$  in automaton  $T_i$  is simulated by  $N \times (d_{max} + 2)$  transitions in  $\mathcal{N}$  as follows. For each  $(z, j) \in \{0, 1, \dots, d_{max}, \infty\} \times \{1, \dots, N\}$ , a transition removes a token from the place  $p_{\ell_i^0}$  corresponding to the unique initial location  $\ell_i^0$ , a token from  $p_{(a,z,j)}^i$  and adds a token to the place  $p_{\ell'}$  corresponding to  $\ell'$ . The deadline is checked on the arc from the place  $p_{(a,z,j)}^i$  by a constraint which checks the age of the token to be in the interval  $[0, z]$ . As any deadline value is possible, and any of the  $N$  places can be filled, one of the  $N \times (d_{max} + 2)$  transitions is non-deterministically chosen.

A transition of the form  $(\ell, g, i!j(a[d]), Y, \ell')$  in automaton  $T_i$  is simulated in a similar way by  $N + 1$  transitions. In each of the  $N$  of these transitions, tokens for control locations are added and removed as in the previous case. For each  $k \in \{1, \dots, N\}$ , one of the  $N$  transitions adds a token to the place  $p_{a,d,k}^j$  if it is empty. The  $(N + 1)$ -th transition simulates the possibility that the task  $a$  is not relevant (only  $N$  are relevant at any point) and so it simulates only the change in control location and adds no other tokens. One of these  $N + 1$  transitions is chosen non-deterministically. Observe that the first  $N$  transitions add a token only to an empty place  $p_{a,d,k}^j$  by definition of an 1-safe Petri net.

Clock resets are simulated by adding and removing a token from the corresponding place  $p_x$  for the clock. Clock constraints are simulated by read arcs. These arcs are connected with the corresponding transitions that are described above.

The formal construction is in [1]. Thus, the control state reachability in  $\mathcal{M}$  to reach  $(f_1, \dots, f_N) \in L_1 \times \dots \times L_N$  reduces to the coverability problem of the marking  $M$  given by  $M(p_{f_i}) = 1$  for all  $1 \leq i \leq N$  (and hence  $M(p_\ell) = 0$  for all  $\ell \notin \{f_1, \dots, f_N\}$ ). The control state reachability of  $\mathcal{M}$  thus reduces to the coverability of the constructed 1-safe timed Petri net with read arcs. Since the coverability of 1-safe timed Petri nets with read arcs is PSPACE-complete [17], the control state reachability of  $\mathcal{M}$  is also PSPACE-complete. ◀

## 5 Configuration Reachability

In this section, we explore the general question of the configuration reachability problem for  $N$ -MTA. We first show (theorem 4) that the configuration reachability problem for  $N$ -MTA is undecidable.

► **Theorem 4.** *The configuration reachability problem for  $N$ -MTA is undecidable. This undecidability holds even in the case of time-independent  $N$ -MTA.*

**Proof.** The proof is done by a reduction from the reachability problem for a 2-counter machine (which is known to be undecidable [15]). The main idea is to construct an 1-MTA whose set of states contains the states of the two counter machine plus some auxiliary states

that are used to simulate the zero tests as we will see later on. The 1-MTA has two types of tasks  $a$  and  $b$ . The number of pending tasks of type  $a$  (resp.  $b$ ) corresponds to the value of the counter  $c_1$  (resp.  $c_2$ ). Furthermore, the 1-MTA has one clock  $x$  that is used to check that no time elapsed when simulating some transitions of the two counter machine.

To simulate an increment transition of the form  $(q, c_1++, q')$  (resp.  $(q, c_2++, q')$ ) of the two counter machine, the 1-MTA proceeds as follows: first it checks that the value of the clock  $x = 0$ , then it will change its state from  $q$  to  $q'$  and finally adds a pending task of type  $a$  (resp.  $b$ ) with zero as its deadline. Observe that we need only one transition to perform all these steps of the simulation of an increment operation.

To simulate a decrement transition of the form  $(q, c_1--, q')$  (resp.  $(q, c_2--, q')$ ) of the two counter machine, the 1-MTA proceeds as follows: first it checks that the value of the clock  $x = 0$ , then it will change its state from  $q$  to  $q'$  and finally consumes a pending task of type  $a$  (resp.  $b$ ). Observe that we need only one transition to perform all these steps of the simulation of an increment operation.

To simulate a zero test transition of the form  $(q, c_1 == 0, q')$  (resp.  $(q, c_2 == 0, q')$ ) of the two counter machine, the 1-MTA proceeds as follows: (i) it checks that the value of the clock  $x = 0$ , (ii) it enters to a loop where it consumes a task of type  $b$  (resp.  $a$ ) and creates a task of the same type but its deadline is now set to one time unit, (iii) it will change its state from  $q$  to  $q'$ , (iv) it checks that the value of the clock  $x$  is still zero, (v) it checks that one time unit has elapsed (ie., checking whether  $x \in [1, 1]$ ) and resets the clock  $x$ , (vi) it enters to a loop where it consumes a task of type  $b$  (resp.  $a$ ) and creates a task of the same type but its deadline is now set to zero, and (vii) it checks that the value of  $x = 0$ . Here the auxiliary states are needed in the simulation of these steps.

Observe that if the 1-MTA reaches the final state with empty set of pending tasks, then all the simulation of the zero tests are performed correctly. Finally, note that the constructed 1-MTA is time independent. ◀

We now focus on the class of stateless and time-independent  $N$ -MTA.

► **Theorem 5.** *The configuration reachability problem for stateless and time-independent  $N$ -MTA is decidable.*

We begin by setting up some notations for the proof.

### Well-quasi-orders and Higman's Lemma

Given a set  $\mathcal{Q}$ , a *quasi-order* on  $\mathcal{Q}$  is a reflexive and transitive relation  $\preceq \subseteq \mathcal{Q} \times \mathcal{Q}$ . An infinite sequence  $(q_1, q_2, \dots)$  in  $\mathcal{Q}$  is said to be *saturating* if there exists indices  $i < j$  such that  $q_i \preceq q_j$ . A quasi-order  $\preceq$  is a *well-quasi-order* (wqo) [12] on  $\mathcal{Q}$  if every infinite sequence in  $\mathcal{Q}$  is saturating. Let  $\sqsubseteq$  be a quasi-order on  $\mathcal{Q}$ . The *induced monotone domination order*  $\preceq$  on  $\mathcal{Q}^*$ , (i.e., the set of finite words over  $\mathcal{Q}$ ) is defined as follows:  $a_1 a_2 \dots a_m \preceq b_1 b_2 \dots b_n$  if there exists a strictly increasing function  $g : \{1, 2, \dots, m\} \rightarrow \{1, 2, \dots, n\}$  such that, for all  $1 \leq i \leq m$ ,  $a_i \sqsubseteq b_{g(i)}$ . It is well-known by *Higman's Lemma* [10] that if  $\sqsubseteq$  is a wqo on  $\mathcal{Q}$ , then the induced domination order  $\preceq$  is also a wqo on  $\mathcal{Q}^*$ . As an example, let  $\Sigma = \{1, 2, \dots, 12\}$  and let  $\mathcal{Q}$  be the power set of  $\Sigma$ . Define  $\sqsubseteq$  on  $\mathcal{Q}$  to be the set inclusion relation.  $\sqsubseteq$  is clearly a wqo since  $\mathcal{Q}$  is finite. The induced monotone domination order  $\preceq$  on  $\mathcal{Q}^*$  is the subword order: for example,  $\{1, 2\}\{3\}\{5, 6, 7\} \preceq \{1, 2, 9\}\{1\}\{3, 11\}\{12\}\{4, 5, 6, 7\}$ .

### Encoding Configurations

We have seen in section 3 that a configuration of an  $N$ -MTA  $\mathcal{M}$  is a tuple  $(\bar{q}, \bar{m})$  where  $\bar{q}$  is the sequence of states in each  $T_i$ ,  $1 \leq i \leq N$ , and  $\bar{m}$  is the tuple of multisets  $(M_1, \dots, M_N)$  corresponding to each  $T_i$ . Given  $(\ell_1, \dots, \ell_N) \in L_1 \times \dots \times L_N$ , we are interested in finding whether the configuration  $\mathbf{c}_{goal} = (\bar{q}, \bar{m})$  is reachable, where  $\bar{q} = (q_1, \dots, q_N)$ ,  $q_i = (\ell_i, \mathbf{0})$  and  $\bar{m} = (\emptyset, \dots, \emptyset)$ . A configuration  $\mathbf{c}$  is called *good* if  $\mathbf{c}_{goal}$  is reachable from  $\mathbf{c}$ . A configuration is *bad* if it is not good. Clearly,  $\mathbf{c}_{goal}$  is reachable in  $\mathcal{M}$  iff some initial configuration  $\mathbf{c}_0$  is good.

We now construct an equivalence relation on  $\mathcal{M}$  by encoding the configurations of  $\mathcal{M}$  as words over a certain alphabet. This will enable us to define a wqo on the resulting transition system. Let  $K$  be the maximal constant used in the clock constraints and deadlines in  $\mathcal{M}$ . Let  $[K] = \{0, 1, \dots, K, \infty\}$ . Let  $\mathbf{reg} = \{r_0, r_1, \dots, r_{2K}\}$  be a finite set of *regions*, where for  $0 \leq i \leq K$ ,  $r_{2i}$  is defined as the singleton  $\{i\}$ , while  $r_{2i+1}$  is defined as the interval  $(i, i+1)$  for  $0 \leq i \leq K-1$ . We also define the region  $r_{2K+1}$  as  $(K, \infty)$ . Let  $\Gamma_1$  be the set  $\mathcal{X} \times \mathbf{reg}$ , and let  $\Gamma_2$  be a multiset over  $\{(a, r, j)_i \mid a \in \Sigma, r \in \mathbf{reg}, j \in [K], 1 \leq i \leq N\}$ . Let  $\Gamma_3$  be the set  $\mathcal{X} \times r_{2K+1}$ , and let  $\Gamma_4$  be a multiset over  $\{(a, r_{2K+1}, j)_i \mid a \in \Sigma, 1 \leq i \leq N, j \in [K]\}$ .

Let  $\Upsilon, \Delta$  respectively be the power sets of  $\Gamma_1 \cup \Gamma_2$  and  $\Gamma_3 \cup \Gamma_4$ . Let  $\mathcal{L} = L_1 \times \dots \times L_N$ . We consider words of the form  $\alpha w(P + \epsilon)$  where  $\alpha \in \mathcal{L}$ ,  $w \in \Upsilon^*$  and  $P \in \Delta$ . Since  $\Upsilon, \Delta, \mathcal{L}$  are finite, they are all clearly well-quasi-ordered by set inclusion, and the set of words of the form  $\alpha w(P + \epsilon)$  is well-quasi-ordered by the induced monotone domination order  $\preceq$ :  $\alpha_1 \rho_1 \dots \rho_m P_1 \preceq \alpha_2 \gamma_1 \dots \gamma_n P_2$  if  $\alpha_1 = \alpha_2$ ,  $P_1 \subseteq P_2$ , and there exists a strictly increasing function  $g: \{1, 2, \dots, m\} \rightarrow \{1, 2, \dots, n\}$  such that for all  $1 \leq i \leq m$ ,  $\rho_i \subseteq \gamma_{g(i)}$ .

We next associate to any configuration  $\mathbf{c}$  of  $\mathcal{M}$ , a canonical word  $W(\mathbf{c}) \in \mathcal{L} \cdot \Upsilon^* \cdot (\Delta + \epsilon)$ . Let  $y_{i,1}, \dots, y_{i,|\mathcal{X}_i|}$  be the set of clocks in  $T_i$ . Given a configuration  $\mathbf{c} = (\bar{q}, \bar{m})$  with  $\bar{q} = ((\ell_1, \nu_1), \dots, (\ell_N, \nu_N))$  and  $\bar{m} = (M_1, \dots, M_N)$ ,  $\bar{q}$  is completely specified by describing for each  $1 \leq i \leq N$ , (i) the locations  $\ell_i$ , (ii) the tuples  $(\alpha_{i,j}, \text{frac}(y_{i,j}))$  (resp.  $\alpha_{ij}$ ) if  $\alpha_{ij} = ((y_{i,j}, \mathbf{reg}(\nu(y_{i,j})))$  is in  $\Gamma_1$  (resp.  $\Gamma_3$ ) and  $1 \leq j \leq |\mathcal{X}_i|$ . Observe that here we use  $\text{frac}(y_{i,j})$  (resp.  $\mathbf{reg}(\nu(y_{i,j}))$ ) to denote the fractional part (resp. the corresponding region) of  $\nu(y_{i,j})$ . The former case keeps track of clocks, their regions as well as the fractional parts of the clock valuations, while in the latter, the value of clock  $y_{i,j}$  is more than  $K$ , (iii) the multi set consisting of tuples  $(\beta_i, \text{frac}(\text{age}(a)))$  (resp.  $\beta_i$ ) if  $\beta_i = (a, \mathbf{reg}(\text{age}(a)), d)$  is in  $\Gamma_2$  (resp.  $\Gamma_4$ ). The former keeps track of tasks, the region of their ages, and their deadlines, along with the fractional parts of the ages, while in the latter, the age of the task is more than  $K$ . Observe that  $\text{age}(a)$  returns the age of the task  $a$ .

Next, we group together the symbols  $\alpha_h \in \Gamma_1, \beta_g \in \Gamma_2$  having the same fractional parts. Notice that the fractional parts are retained only for clocks (tasks) whose value (age) has not yet exceeded  $K$ . This yields a new set of  $\Gamma_1 \cup \Gamma_2$  letters paired with their fractional parts  $\{(\zeta_i, \text{frac}_i) \mid 1 \leq i \leq p\}$  where  $\zeta_i$  is a (multi)set of symbols from  $\Gamma_1 \cup \Gamma_2$  and  $\text{frac}_i$  is the fractional part of those symbols.  $p$  is the number of distinct fractional parts in  $\mathbf{c}$ . We then form the word  $w = \rho_{i_{z_1}} \dots \rho_{i_{z_p}} \in \Upsilon^+$  where  $z_1 \dots z_p$  is a permutation of  $1 \dots p$  that puts  $\text{frac}_{z_1} \dots \text{frac}_{z_p}$  in ascending order. Let  $P \in \Delta$  be the set obtained (if any) by grouping all the symbols  $\alpha_h \in \Gamma_3$  and  $\beta_g \in \Gamma_4$ . We then define  $W(\mathbf{c}) = \alpha.w.P \in \mathcal{L} \cdot \Upsilon^* \cdot (\Delta + \epsilon)$  as the canonical word encoding  $\mathbf{c}$ .

► **Example 6.** Consider a 2-MTA  $\mathcal{M}$ . Let  $x_1, x_2$  be the clocks of  $T_1$  and  $y_1, y_2$  be the clocks of  $T_2$ . Let  $K = 3$  be the maximal constant used in  $\mathcal{M}$ . Consider the configurations  $\mathbf{c}_1 = ((s_1, 0.5, 2.1), (s_2, 1.7, 2.5), (\{(a, 1.1, 2), (b, 2.3, \infty), (c, 3.5, \infty)\}, \{(d, 1.9, 2), (e, 0.7, 1)\}))$  and  $\mathbf{c}_2 = ((s_1, 0.5, 2.4), (s_2, 1.9, 2.5), (\{(a, 1.4, 2), (b, 2.45, \infty), (c, 3.9, \infty)\}, \{(d, 1.99, 2), (e, 0.9, 1)\}))$ .

Then  $W(c_1) = W(c_2) = \alpha w P$  where  $\alpha = (s_1, s_2)$ ,  $P = \{(c, r_7, \infty)_1\}$ , and  $w = \{(x_2, r_5), (a, r_3, 2)_1\} \{(b, r_5, \infty)_1\} \{(x_1, r_1), (y_2, r_5)\} \{(y_1, r_3), (e, r_1, 1)_2\} \{(d, r_3, 2)_2\}$ .

Two configurations  $c_1, c_2$  are equivalent ( $c \sim c'$ ) if  $W(c_1) = W(c_2)$ . A configuration  $c_1$  is dominated by a configuration  $c_2$  (written  $c_1 \preceq c_2$ ) if writing  $c_2 = (\bar{q}_2, \bar{m}_2)$ , there exists  $\bar{q}_1, \bar{m}_1$  such that  $\bar{m}_1 = (M'_1, \dots, M'_N)$  with  $M'_i \subseteq M_i$  for all  $i$ , and  $c_1 \sim (\bar{q}_1, \bar{m}_1)$ . It can be easily seen that  $c_1 \preceq c_2$  iff  $W(c_1) \preceq W(c_2)$ . In fact, the following lemma shows that  $\sim$  is a bisimulation relation.

► **Lemma 7.** *Let  $c_1, c_2$  be two configurations of an  $N$ -MTA. Let  $e \in E_i$  be a transition,  $1 \leq i \leq N$ , and let  $t \in \mathbb{R}$ . If  $c_1 \sim c_2$ , then*

- (1) *If  $c_1 \xrightarrow{e} c'_1$ , there exists  $c'_2$  such that  $c_2 \xrightarrow{e} c'_2$  and  $c'_1 \sim c'_2$ . If  $c_2 \xrightarrow{e} c'_2$ , there exists  $c'_1$  such that  $c_1 \xrightarrow{e} c'_1$  and  $c'_1 \sim c'_2$ .*
- (2) *If  $c_1 \xrightarrow{t} c'_1$ , there exists  $c'_2$  and  $t' \in \mathbb{R}$  such that  $c_2 \xrightarrow{t'} c'_2$  and  $c'_1 \sim c'_2$ . If  $c_2 \xrightarrow{t} c'_2$ , there exists  $c'_1$  and  $t' \in \mathbb{R}$  such that  $c_1 \xrightarrow{t'} c'_1$  and  $c'_1 \sim c'_2$ .*

As an easy corollary of the above, we see that  $\sim$  preserves goodness and badness: For any configurations  $c \sim c'$ ,  $c$  is good iff  $c'$  is good. The proof follows from the definition of goodness and Lemma 7, whose proof can be found in the extended version of the paper [1].

It is hence sufficient to only consider configurations upto  $\sim$ -equivalence, and we define the quotient labeled transition system  $\mathcal{M}/\sim$  to consist of all the words  $W(c)$  whenever  $c$  is a configuration of  $\mathcal{M}$ . Call  $\mathcal{M}/\sim$  as  $\mathcal{W}$ .  $\mathcal{W} = \{W(c) \mid c \text{ is a configuration in } \mathcal{M}\}$ . For  $W_1, W_2 \in \mathcal{W}$ , and a transition  $e \in E_i$  for  $1 \leq i \leq N$ , we define a transition  $W_1 \xrightarrow{e} W_2$  if for all  $c_1 \in W^{-1}(W_1)$ , there is some configuration  $c_2 \in W^{-1}(W_2)$  such that  $c_1 \xrightarrow{e} c_2$ . The timed transition is defined similarly. Corresponding to each initial configuration  $c_0$  in  $\mathcal{M}$ , we consider  $W_0 = W(c_0)$  to be an initial word in  $\mathcal{W}$ . Let  $\mathcal{W}_0$  be the set of initial words corresponding to initial configurations  $c_0$ . It can be seen that for any  $W_1, W_2 \in \mathcal{W}$ , and a transition  $e \in E_i$  or  $t \in \mathbb{R}$ ,  $W_1 \xrightarrow{\alpha} W_2$  ( $\alpha \in \{e, t\}$ ) iff there exist configurations  $c_1 \in W^{-1}(W_1)$  and  $c_2 \in W^{-1}(W_2)$  such that  $c_1 \xrightarrow{\alpha} c_2$ . Given a word  $W \in \mathcal{W}$ , and  $\alpha \in \{e, t\}$  for some transition  $e$  and time  $t \in \mathbb{R}$ , let  $\text{succ}(W) = \{W' \in \mathcal{W} \mid W \xrightarrow{\alpha} W'\}$  denote the successors of  $W$  in  $\mathcal{W}$ .

► **Lemma 8.** *For any word  $W$ , the set  $\text{succ}(W)$  is finite and effectively computable.*

Let  $\mathcal{W}_0$  be the set of initial words corresponding to  $W(c_0)$  for initial configurations  $c_0$ . Let  $W_\emptyset = W(c_{\text{goal}})$ . Algorithm 1 decides whether the configuration  $c_{\text{goal}}$  can be reached. In this algorithm, the function  $\text{Minimize}(R)$  is used, where  $R \subseteq \mathcal{W}$  is a set of words. It does the following: it chooses a word  $W_1 \in R$  and removes  $W_1$  from  $R$  if there exists a word  $W_2 \in R$  such that  $W_2 \preceq W_1$ , and then repeats the procedure until all words in  $R$  are processed. Overall, the algorithm works as follows. Till the set **Next** of words waiting to be processed is non-empty, the algorithm chooses one word from **Next**, and moves it to the **Processed** set. It also generates all successors of the chosen word, minimizes them, and adds them to **Next** unless there is already some  $\preceq$ -smaller word in **Next** or **Processed**. If a new word is added to **Next**, the algorithm removes at the same time all  $\preceq$ -bigger words from both **Next** and **Processed**. The correctness of the algorithm is discussed next.

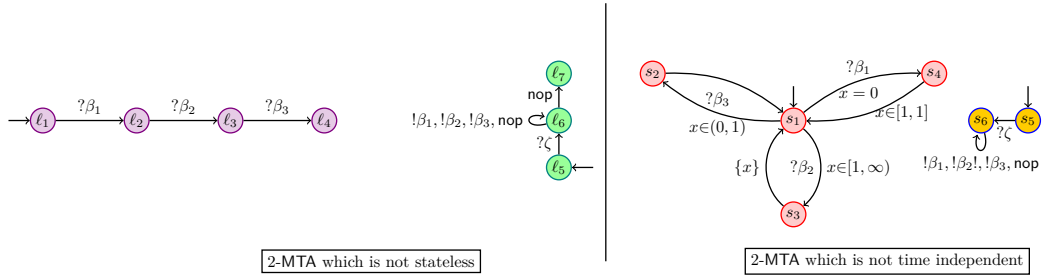
A set of words  $R$  is good (denoted  $\text{Good}(R)$ ) iff there exists some word  $W \in R$  which is good. A word  $W$  is good iff there exists a good configuration  $c$  such that  $W(c) = W$ . If  $W$  is a good word, and if  $i \in \mathbb{N}$  is the length of the shortest path (excluding time elapse transitions) from  $W$  to  $W_\emptyset$ , then we say that  $\text{dist}(W)$  is  $i$ . Given a set  $R$  of words,  $\text{dist}(R) \in \mathbb{N} \cup \{\infty\}$  is defined as the length of the shortest path (excluding time elapse transitions) from some  $W \in R$  to  $W_\emptyset$ . More precisely, if  $R = \emptyset$ , then  $\text{dist}(R) = \infty$ , otherwise,  $\text{dist}(R) = \min_{W \in R} \text{dist}(W)$ .

**Algorithm 1** REACH EMPTY.

---

**Input:** A stateless, time-independent  $N$ -MTA, and configuration  $\mathbf{c}_{goal} = (\bar{q}, \bar{m})$  as above.  
**Output:** TRUE if  $\mathbf{c}_{goal}$  is reachable. Otherwise, FALSE.  
**if**  $W_\emptyset \in \mathcal{W}_0$ , **then return** TRUE;  
 Processed =  $\emptyset$ ;  
 Next = Minimize( $\mathcal{W}_0$ );  
**while** Next  $\neq \emptyset$  **do**  
   **leftmargin=0.5in** Pick and remove a word  $W$  from Next and move it to Processed,  
   **leftmargin=0.5iin** **foreach**  $U \in \text{Minimize}(\text{succ}(W))$   
     **leftmargin=0.5in** **if**  $U = W_\emptyset$ , **then return** TRUE,  
     **leftmargin=0.5iin** **else if**  $\nexists V \in \text{Processed} \cup \text{Next}$  s.t.  $V \preceq U$ ,  
       **then**  
     **leftmargin=0.5iin** Remove all  $V$  from  $\text{Processed} \cup \text{Next}$  s.t.  $U \preceq V$   
     **leftmargin=0.5ivn** Add  $U$  to Next  
**return** FALSE

---

■ **Figure 3**► **Lemma 9.**

1.  $\text{Good}(\text{Processed} \cup \text{Next}) \rightarrow \text{Good}(\mathcal{W}_0)$
2.  $\text{Good}(\mathcal{W}_0) \rightarrow \text{dist}(\text{Processed}) > \text{dist}(\text{Next})$

To prove the invariants, we use the following lemma.

► **Lemma 10.** *If  $W \preceq W'$  and  $\text{dist}(W') = i$ , then  $\text{dist}(W) = j$  for some  $j \leq i$ .*

Due to the well-quasi ordering, the algorithm terminates: if not, over a period of time, there will be an infinite sequence of words in Next, each new word added having the property that it does not dominate any of its predecessors. This would constitute an infinite non saturating sequence, directly contradicting Higman's Lemma. The algorithm returns FALSE only when Next is empty. Then,  $\text{dist}(\text{Processed}) > \text{dist}(\text{Next})$  is not true. Therefore, by invariant 2 in lemma 9,  $\mathcal{W}_0$  is not good. The algorithm returns TRUE only if either  $W_\emptyset$  is already in  $\mathcal{W}_0$ , or if  $W_\emptyset$  is a member of  $\text{Minimize}(\text{succ}(W))$  for some  $W \in \text{Next}$ . In either case, Next is good. Then, by invariant 1 of lemma 9,  $\mathcal{W}_0$  is good. This gives the following lemma.

► **Lemma 11.** *Algorithm REACH EMPTY terminates and returns true iff starting from the initial configuration  $\mathbf{c}_0$  in  $\mathcal{M}$ ,  $\mathbf{c}_{goal}$  is reachable.*

This concludes the proof of theorem 5. A detailed discussion with proofs for the lemmas can be found in the extended version of the paper [1].

Notice that the stateless, and time-independent properties of  $\mathcal{M}$  are crucial in Lemma 10. The example below shows that relaxing either condition violates Lemma 10.



To the left is a 2-MTA which is not stateless. It can be seen that  $c_1 = (\ell_1, \ell_6, \{\beta_1, \beta_3\}, \emptyset) \preceq (\ell_1, \ell_6, \{\beta_1, \beta_2, \beta_3\}, \emptyset) = c_2$ . Hence,  $W(c_1) \preceq W(c_2)$ . Indeed from  $c_2$ , one can reach  $(\ell_4, \ell_6, \emptyset, \emptyset)$ , but not from  $c_1$ . To the right is a 2-MTA which is not time independent. It can be seen that  $c_1 = (((s_1, 0), s_6), \{(\beta_1, 0, \infty), (\beta_3, 0, \infty)\}, \emptyset) \preceq (((s_1, 0), s_6), \{(\beta_1, 0, \infty), (\beta_2, 0, \infty), (\beta_3, 0, \infty)\}, \emptyset) = c_2$ . However,  $((s_1, 0), s_6), \emptyset, \emptyset$  is reachable from  $c_2$  but not from  $c_1$ .

## 6 Conclusion

We proposed a model to address the verification problem for timed asynchronous programs. We identified a special subclass (stateless and time-independent) for which the reachability problem is decidable and control reachability is PSPACE-complete. There are multiple avenues for further work. The first question is to check the tightness of the EXPSPACE lower bound provided. Another question would be to consider the model where we use *priority bags* instead of bags. In a priority bag, tasks have associated deadlines and priorities. The process, while picking up a task for execution, is expected to pick up a task with the highest priority. Queues are yet another interesting data structure in place of bags: in this set up, the tasks which require a processor's attention are picked up in the order in which they were assigned by various processes. One can also look at mutiset timed pushdown systems, which extend the model of [16] with time, and multiple processes. Finally, we can move from the one player setting to two players, where the environment chooses a task for the process to execute. Under this two player setting, the question would be if the system has a strategy to execute all the pending tasks.

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